

Foundation of Cryptography (0368-4162-01), Lecture 1

One Way Functions

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Section 1

Notation

Notation I

- For $t \in \mathbb{N}$, let $[t] := \{1, \dots, t\}$.
- Given a string $x \in \{0, 1\}^*$ and $0 \leq i < j \leq |x|$, let $x_{i, \dots, j}$ stands for the substring induced by taking the i, \dots, j bit of x (i.e., $x[i] \dots, x[j]$).
- Given a function f defined over a set \mathcal{U} , and a set $\mathcal{S} \subseteq \mathcal{U}$, let $f(\mathcal{S}) := \{f(x) : x \in \mathcal{S}\}$, and for $y \in f(\mathcal{U})$ let $f^{-1}(y) := \{x \in \mathcal{U} : f(x) = y\}$.
- poly stands for the set of all polynomials.
- The worst-case running-time of a *polynomial-time algorithm* on input x , is bounded by $p(|x|)$ for some $p \in \text{poly}$.
- A function is *polynomial-time computable*, if there exists a polynomial-time algorithm to compute it.

Notation II

- PPT stands for probabilistic polynomial-time algorithms.
- A function $\mu: \mathbb{N} \mapsto [0, 1]$ is negligible, denoted $\mu(n) = \text{neg}(n)$, if for any $p \in \text{poly}$ there exists $n' \in \mathbb{N}$ with $\mu(n) \leq 1/p(n)$ for any $n > n'$.

Distribution and random variables I

- The support of a distribution P over a finite set \mathcal{U} , denoted $\text{Supp}(P)$, is defined as $\{u \in \mathcal{U} : P(u) > 0\}$.
- Given a distribution P and an event E with $\Pr_P[E] > 0$, we let $(P \mid E)$ denote the conditional distribution P given E (i.e., $(P \mid E)(x) = \frac{P(x) \wedge E}{\Pr_P[E]}$).
- For $t \in \mathbb{N}$, let U_t denote a random variable uniformly distributed over $\{0, 1\}^t$.
- Given a random variable X , we let $x \leftarrow X$ denote that x is distributed according to X (e.g., $\Pr_{x \leftarrow X}[x = 7]$).
- Given a finite set S , we let $x \leftarrow S$ denote that x is uniformly distributed in S .

Distribution and random variables II

- We use the convention that when a random variable appears twice in the same expression, it refers to a *single* instance of this random variable. For instance, $\Pr[X = X] = 1$ (regardless of the definition of X).
- Given distribution P over \mathcal{U} and $t \in \mathbb{N}$, we let P^t over \mathcal{U}^t be defined by $D^t(x_1, \dots, x_t) = \prod_{i \in [t]} D(x_i)$.
- Similarly, given a random variable X , we let X^t denote the random variable induced by t independent samples from X .

Section 2

One Way Functions

One-Way Functions

Definition 1 (One-Way Functions (OWFs))

A polynomial-time computable function $f : \{0, 1\}^* \mapsto \{0, 1\}^*$ is one-way, if for any PPT A

$$\Pr_{y \leftarrow f(U_n)}[A(1^n, y) \in f^{-1}(y)] = \text{neg}(n)$$

U_n : a random variable uniformly distributed over $\{0, 1\}^n$

polynomial-time computable: there exists a polynomial-time algorithm F , such that $F(x) = f(x)$ for every $x \in \{0, 1\}^*$

PPT: probabilistic polynomial-time algorithm

neg: a function $\mu: \mathbb{N} \mapsto [0, 1]$ is a *negligible* function of n , denoted $\mu(n) = \text{neg}(n)$, if for any $p \in \text{poly}$ there exists $n' \in \mathbb{N}$ such that $\mu(n) < 1/p(n)$ for all $n > n'$

We will typically omit 1^n from the parameter list of A

- ❶ Is this the right definition?
 - Asymptotic
 - Efficiently computable
 - On the average
 - Only against PPT's
- ❷ (most) Crypto implies OWFs
- ❸ Do OWFs imply Crypto?
- ❹ Where do we find them
- ❺ Non uniform OWFs

Definition 2 (Non-uniform OWF))

A polynomial-time computable function $f : \{0, 1\}^* \mapsto \{0, 1\}^*$ is one-way, if for any polynomial-size family of circuits $\{C_n\}_{n \in \mathbb{N}}$

$$\Pr_{y \leftarrow f(U_n)}[C_n(y) \in f^{-1}(y)] = \text{neg}(n)$$

Length preserving functions

Definition 3 (length preserving functions)

A function $f : \{0, 1\}^* \mapsto \{0, 1\}^*$ is length preserving, if $|f(x)| = |x|$ for any $x \in \{0, 1\}^*$

Theorem 4

Assume that OWFs exist, then there exist length-preserving OWFs

Proof idea: use the assumed OWF to create a length preserving one

Partial domain functions

Definition 5 (Partial domain functions)

For $m, \ell: \mathbb{N} \mapsto \mathbb{N}$, let $h: \{0, 1\}^{m(n)} \mapsto \{0, 1\}^{\ell(n)}$ denote a function defined over input lengths in $\{m(n)\}_{n \in \mathbb{N}}$, and maps strings of length $m(n)$ to strings of length $\ell(n)$.

The definition of one-wayness naturally extends to such functions.

OWFs imply Length Preserving OWFs cont.

Let f be a OWF, let $p \in \text{poly}$ be a bound on its computing-time and assume wlg. that p is monotonly increasing (can we?).

Construction 6 (the length preserving function)

Define $g: \{0, 1\}^{p(n)} \mapsto \{0, 1\}^{p(n)}$ as

$$g(x) = f(x_1, \dots, x_n), 0^{p(n) - |f(x_1, \dots, x_n)|}$$

Note that g is length preserving and efficient (why?).

Claim 7

g is one-way.

How can we prove that g is one-way?

Answer: using reduction

Proving that g is one-way

Proof:

Assume that g is not one-way. Namely, there exists PPT A a $q \in \text{poly}$ and an infinite $\mathcal{I} \subseteq \{p(n) : n \in \mathbb{N}\}$, with

$$\Pr_{y \leftarrow g(U_n)}[A(y) \in g^{-1}(y)] > 1/q(n) \quad (1)$$

for any $n \in \mathcal{I}$ (?).

We would like to use A for inverting f .

Algorithm 8 (The inverter B)

Input: 1^n and $y \in \{0, 1\}^*$.

- 1 Let $x = A(1^{p(n)}, y, 0^{p(n)-|y|})$.
- 2 Return $x_{1,\dots,n}$.

Claim 9

Let $\mathcal{I}' := \{n \in \mathbb{N} : p(n) \in \mathcal{I}\}$. Then

- 1 \mathcal{I}' is infinite
- 2 For any $n \in \mathcal{I}'$, it holds that $\Pr_{y \leftarrow g(U_n)}[B(y) \in f^{-1}(y)] > 1/q(p(n))$.

in contradiction to the assumed one-wayness of f . \square

Conclusion

Remark 10

- We directly related the hardness of f to that of g
- The reduction is not “security preserving”

From partial domain functions to all-length functions

Construction 11

Given a function $f: \{0, 1\}^{m(n)} \mapsto \{0, 1\}^{\ell(n)}$,
 $f_{all}: \{0, 1\}^* \mapsto \{0, 1\}^*$ as

$$f_{all}(x) = f(x_1, \dots, x_{k(n)}), 0^{n-k(n)}$$

where $n = |x|$ and $k(n) := \max\{m(n') \leq n: n' \in \mathbb{N}\}$.

Claim 12

Assume that f is a one-way function and that m is monotone, polynomial-time computable and satisfies $\frac{m(n+1)}{m(n)} \leq p(n)$ for some $p \in \text{poly}$, then f_{all} is a one-way function. Further, if f is length preserving, then so is f_{all} .

Proof: ?

Weak One Way Functions

Definition 13 (weak one-way functions)

A polynomial-time computable function $f : \{0, 1\}^* \mapsto \{0, 1\}^*$ is α -one-way, if

$$\Pr_{y \leftarrow f(U_n)}[A(1^n, y) \in f^{-1}(y)] \leq \alpha(n)$$

for any PPT A and large enough $n \in \mathbb{N}$.

- 1 (strong) OWF according to Definition 1, are $\text{neg}(n)$ -one-way according to the above definition
- 2 Examples
- 3 Can we “amplify” weak OWF to strong ones?

Strong to weak OWFs

Claim 14

Assume there exists OWFs, then there exist functions that are $\frac{1}{3}$ one-way, but not (strong) one-way

Proof: let f be a owf. Define $g(x) = (1, g(x))$ if $x_1 = 1$, and 0 otherwise.

Weak to Strong OWFs

Theorem 15

Assume there exists $(1 - \alpha)$ -weak OWFs with $\alpha(n) > 1/p(n)$ for some $p \in \text{poly}$, then there exists (strong) one-way functions.

Proof: we assume wlg that f is length preserving (can we do so?)

Construction 16 (g – the strong one-way function)

Let $t: \mathbb{N} \mapsto \mathbb{N}$ be a polynomial-time computable function satisfying $t(n) \in \omega(\log n / \alpha(n))$. Define $g: (\{0, 1\}^n)^{t(n)} \mapsto (\{0, 1\}^n)^{t(n)}$ as

$$g(x_1, \dots, x_t) = f(x_1), \dots, f(x_t)$$

Claim 17

g is one-way.

Proving that g is one-way – the naive approach

Let A be a potential inverter for g , and assume that A tries to attacks each of the t outputs of g *independently*. Then

$$\Pr_{y \leftarrow g(U_n^{t(n)})} [A(y) \in g^{-1}(y)] \leq (1 - \alpha(n))^{t(n)} \leq e^{-\omega(\log n)} = \text{neg}(n)$$

A less naive approach would be to assume that A goes over output sequentially.

Unfortunately, we can assume none of the above.

Failing Sets

Definition 18 (failing set)

A function $f : \{0, 1\}^n \mapsto \{0, 1\}^{\ell(n)}$ has a $(\delta(n), \varepsilon(n))$ -failing set for A , if for large enough n , exists set $\mathcal{S}(n) \subseteq \{0, 1\}^{\ell(n)}$ with

- 1 $\Pr[f(U_n) \in \mathcal{S}(n)] \geq \delta(n)$, and
- 2 $\Pr[A(y) \in f^{-1}(y)] < \varepsilon(n)$, for every $y \in \mathcal{S}(n)$

Claim 19

Let f be a $(1 - \alpha)$ -OWF. Then f has $(\alpha(n)/2, 1/p(n))$ -failing set for any PPT A and $p \in \text{poly}$.

Proof: Assume \exists PPT A , a $p \in \text{poly}$ and an infinite set $\mathcal{I} \subseteq \mathbb{N}$ such that for every $n \in \mathcal{I}$, $\exists \mathcal{L}(n) \subseteq \{0, 1\}^n$ with

- 1 $\Pr[f(U_n) \in \mathcal{L}(n)] \geq 1 - \alpha(n)/2$, and
- 2 $\Pr[A(y) \in f^{-1}(y)] \geq 1/p(n)$, for every $y \in \mathcal{L}(n)$

We'll use A to contradict the hardness of f .

Using A to invert f

Algorithm 20 (The inverter B)

Input: $y \in \{0, 1\}^n$.

Do (with fresh randomness) for $np(n)$ times:

If $x = A(y) \in f^{-1}(y)$, return x

Clearly, B is a PPT

Claim 21

For every $n \in \mathcal{I}$, it holds that

$$\Pr_{y \leftarrow f(U_n)}[B(y) \in f^{-1}(y)] > 1 - \alpha(n)$$

Hence, f is not $(1 - \alpha(n))$ -one-way \square

Proof of Claim 21(all probabilities below are also over $y \leftarrow f(U_n)$):

$$\begin{aligned} & \Pr[B(y) \in f^{-1}(y)] \\ & \geq \Pr[B(y) \in f^{-1}(y) \wedge y \in \mathcal{L}(n)] \\ & = \Pr[y \in \mathcal{L}(n)] \cdot \Pr[B(y) \in f^{-1}(y) \mid y \in \mathcal{L}(n)] \\ & \geq (1 - \alpha(n)/2) \cdot (1 - (1 - 1/p(n))^{np(n)}) \\ & \geq (1 - \alpha(n)/2) \cdot (1 - 2^{-n}) > 1 - \alpha(n). \square \end{aligned}$$

Proving that g is one-way

We show that if g is not OWF, then f has no flailing-set of the “right” type.

Claim 22

Assume \exists PPT A , $p \in \text{poly}$ and an infinite set $\mathcal{I} \subseteq \mathbb{N}$ s.t.

$$\Pr_{z \leftarrow g(U_n^{t(n)})}[A(z) \in g^{-1}(z)] \geq 1/p(n) \quad (2)$$

for every $n \in \mathcal{I}$. Then \exists PPT B and $q \in \text{poly}$ s.t.

$$\Pr_{y \leftarrow \mathcal{S}}[B(y) \in f^{-1}(y)] \geq 1/q(n) \quad (3)$$

for every $n \in \mathcal{I}$ and $\mathcal{S} \subseteq \{0, 1\}^n$ with $\Pr_{y \leftarrow f(U_n)}[\mathcal{S}] \geq \alpha(n)/2$.

Namely, f does not have a $(\alpha(n)/2, 1/q(n))$ -failing set.

Algorithm B

Algorithm 23 (No failing set algorithm B)

Input: $y \in \{0, 1\}^n$.

- 1 Choose $z = (z_1, \dots, z_t) \leftarrow g(U_n^t)$ and $i \leftarrow [t]$
- 2 Set $z' = (z_1, \dots, z_{i-1}, y, z_{i+1}, \dots, z_t)$
- 3 Return $A(z')_i$

Fix $n \in \mathcal{I}$ and a set $\mathcal{S} \subseteq \{0, 1\}^n$ of the right probability. We analyze B's success probability using the following (inefficient) algorithm B*:

Algorithm B*

Definition 24 (Bad)

For $z \in \text{Im}(g)$ (the image of g), we set $\text{Bad}(z) = 1$ iff $\nexists i \in [t]$ with $z_i \in \mathcal{S}$.

B^* differ from B in the way it chooses z' : in case $\text{Bad}(z) = 1$, it sets $z' = z$. Otherwise, it sets i to an arbitrary index $j \in [t]$ with $z_j \in \mathcal{S}$, and sets z' as B does with respect to this i .

Claim 25

$$\Pr_{y \leftarrow \mathcal{S}}[B^*(y) \in f^{-1}(y)] \geq \frac{1/p(n)}{-} \text{neg}(n),$$

and therefore $\Pr_{y \leftarrow \mathcal{S}}[B(y) \in f^{-1}(y)] \geq \frac{1}{t(n)p(n)} - \text{neg}(n). \square$

Claim 25 follows from the following two claims,

Claim 26

$$\Pr_{z \leftarrow g(U_n^t)}[\text{Bad}(z)] = \text{neg}(n)$$

Claim 27

Let $Z = g(U_n^t)$ and let Z' be the value of z' induced by a random execution of B^* on $y \leftarrow (f(U_n) \mid f(U_n) \in S)$. Then Z and Z' are identically distributed.

The claims imply Claim 25.

$$\Pr_{y \leftarrow \mathcal{S}}[B^*(y) \in f^{-1}(y)] \geq \Pr_{z \leftarrow g(U_n^t)}[A(z) \in g^{-1}(z) \wedge \neg \text{Bad}(z)] \quad (4)$$

$$\begin{aligned} & \Pr_{z \leftarrow g(U_n^t)}[A(z) \in g^{-1}(z)] \\ & \leq \Pr[A(z) \in g^{-1}(Z) \wedge \neg \text{Bad}(z)] + \Pr[\text{Bad}(z)] \end{aligned} \quad (5)$$

It follows that

$$\begin{aligned} \Pr_{y \leftarrow \mathcal{S}}[B^*(y) \in f^{-1}(y)] & \geq \Pr_{z \leftarrow g(U_n^t)}[A(z) \in g^{-1}(z)] - \text{neg}(n) \\ & \geq \frac{1}{p(n)} - \text{neg}(n). \square \end{aligned}$$

Proof of Claim 26?

Proof of Claim 27: Consider the following process for sampling Z_i :

- 1 Let $\beta = \Pr_{y \leftarrow f(U_n)}[S]$. Set $\ell_i = 1$ wp β and $\ell_i = 0$ otherwise.
- 2 If $\ell_i = 1$, let $y \leftarrow (f(U_n) \mid f(U_n) \in S)$. Otherwise, set $y \leftarrow (f(U_n) \mid f(U_n) \notin S)$.

It is easy to see that the above process is correct (samples Z correctly).

Now all that B^* does, is repeating Step 2 for one of the i 's with $\ell_i = 1$ (if such exists) \square

Conclusion

Remark 28 (hardness amplification via parallel repetition)

- Can we give a more efficient (secure) reduction?
- Similar theorems for other cryptographic primitives (e.g., Captchas, general protocols)?

What properties of the weak OWF have we used in the proof?